

A CHARACTERIZATION FOR DECIDABLE SEPARABILITY BY PIECEWISE TESTABLE LANGUAGES

WOJCIECH CZERWIŃSKI, WIM MARTENS, LORIJN VAN ROOIJEN, MARC ZEITOUN,
AND GEORG ZETZSCHE

ABSTRACT. The separability problem for word languages of a class \mathcal{C} by languages of a class \mathcal{S} asks, for two given languages I and E from \mathcal{C} , whether there exists a language S from \mathcal{S} that includes I and excludes E , that is, $I \subseteq S$ and $S \cap E = \emptyset$. In this work, we assume some mild closure properties for \mathcal{C} and study for which such classes \mathcal{C} , separability by piecewise testable languages (PTL) is decidable. We characterize these classes in terms of decidability of (two variants of) an unboundedness problem. From this we deduce that separability by PTL is decidable for a number of language classes, such as the context-free languages and languages of labeled vector addition systems. Furthermore, it follows that separability by PTL is decidable if and only if one can compute for any language of the class its downward closure wrt. the scattered substring ordering (i.e., if the set of scattered substrings of any language of the class is effectively regular).

The obtained decidability results contrast some undecidability results. In fact, for all the (non-regular) language classes we present as examples with decidable separability, it is undecidable whether a given language is a PTL itself.

Our characterization involves a result of independent interest, which states that for *any* kind of languages I and E , non-separability is equivalent to the existence of common patterns in I and E .

1. INTRODUCTION

We say that language I can be *separated* from E by language S if S includes I and excludes E , that is, $I \subseteq S$ and $S \cap E = \emptyset$. In this case, we call S a *separator*. We study the *separability problem* of a class \mathcal{C} by a class \mathcal{S} :

Given: Two languages I and E from a class \mathcal{C} .

Question: Can I and E be separated by some language from \mathcal{S} ?

Separability is a classical problem in mathematics and computer science that recently found much new interest. For example, recent work investigated the separability problem of regular languages by piecewise testable languages (Place et al., 2013b; Czerwiński et al., 2013b), locally testable and locally threshold testable languages (Place et al., 2013a) or by first order definable languages (Place and Zeitoun, 2014b). Another recent example is the proof of Leroux (Leroux, 2010) for the decidability of reachability for vector addition systems or Petri nets. Leroux' proof uses separation and goes beyond regular languages. It greatly simplifies earlier proofs by Mayr (Mayr, 1984) and Kosaraju (Kosaraju, 1982).

In this paper we focus on the theoretical underpinnings of separation by piecewise testable languages. Our interest in piecewise testable languages is mainly because of the following two reasons. First, it was shown recently (Place et al., 2013b; Czerwiński et al., 2013b) that separability

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of regular languages (given by their non-deterministic automata) by piecewise testable languages is in PTIME. We found the tractability of this problem to be rather surprising.

Second, piecewise testable languages form a very natural class in the sense that they only reason about the *order* of symbols. More precisely, they are finite Boolean combinations of regular languages of the form $A^*a_1A^*a_2A^*\dots A^*a_nA^*$ in which $a_i \in A$ for every $i = 1, \dots, n$. We are investigating to which extent piecewise testable languages and fragments thereof can be used for computing simple explanations for the behavior of complex systems (Hofman and Martens, 2015).

Separation and Characterization. For classes \mathcal{C} effectively closed under complement, separation of \mathcal{C} by \mathcal{S} is a natural generalization of the *characterization problem* of \mathcal{C} by \mathcal{S} , which is defined as follows. For a given language $L \in \mathcal{C}$ decide whether L is in \mathcal{S} . Indeed, L is in \mathcal{S} if and only if L can be separated from its complement by a language from \mathcal{S} . The characterization problem is well studied. The starting points were famous works of Schützenberger (Schützenberger, 1965) and Simon (Simon, 1975), who solved it for the regular languages by the first-order definable languages and piecewise testable languages, respectively. There were many more results showing that, for regular languages and subclasses thereof (often characterized by a given logic), the problem is decidable, see for example (Brzozowski and Simon, 1971; Zalcstein, 1972; McNaughton, 1974; Knast, 1983; Arfi, 1987; Pin and Weil, 1997; Straubing, 1988; Thérien and Wilke, 1998; Glaßer and Schmitz, 2000; Tesson and Thérien, 2002; Klíma, 2011; Czerwiński et al., 2013a; Place and Zeitoun, 2014a; Place, 2015b; Almeida et al., 2015). Similar problems have been considered for trees (Bojańczyk and Idziaszek, 2009; Benedikt and Segoufin, 2009; Bojańczyk et al., 2012; Antonopoulos et al., 2012; Place and Segoufin, 2015).

Obtaining a decidable characterization for a class is considered as a way to get a fine understanding of the class. In particular, solving a characterization problem requires a proof that a language satisfying some decidable property belongs to the class, which usually yields a canonical construction for the language. For instance, in the case of a logical class, this gives a canonical sentence defining any input language that fulfills the condition we want to prove as a decidable characterization. Recently, several generalizations of the characterization problem have been introduced and investigated. Such generalizations amount to approximating languages of \mathcal{C} by languages of \mathcal{S} . They have been introduced as a mean to obtain even more information on the class \mathcal{S} under study, to then lift this understanding for tackling more complex classes built on top of \mathcal{S} . Separation is the simplest of such approximation problems.

Beyond regular languages. To the best of our knowledge, all the above work and in general all the decidable characterizations were obtained in cases where \mathcal{C} is the class of regular languages, or a subclass of it. This could be due to several negative results which may seem to form a barrier for any nontrivial decidability beyond regular languages. In this work, we consider language classes beyond the regular languages, but we assume them to be full trios (meaning they satisfy some mild closure properties). Beyond the regular languages, we quickly encounter undecidability of the problems above. For instance, for a context-free language (given by a grammar or a pushdown automaton) it is well known that it is undecidable to determine whether it is a regular language, by Greibach’s theorem Greibach (1968). In the same way, one can show that for every full trio that contains the language $\{a^n b^n \mid n \geq 0\}$, it is undecidable to determine whether a given language of the full trio is piecewise testable (see Section 6).

In the case of context-free languages, there is a strong connection between the intersection emptiness problem and separability. Trivially, testing intersection emptiness of two given context-free languages is the same as deciding if they can be separated by some context-free language. However, in general, the negative result is even more overwhelming. Szymanski and Williams (Szymanski and Williams, 1976) proved that separability of context-free languages by regular languages is undecidable. This

was then generalized by Hunt (Hunt III, 1982), who proved that separability of context-free languages by any class containing all the *definite languages* is undecidable. A language L is *definite* if it can be written as $L = F_1 A^* \cup F_2$, where F_1 and F_2 are finite languages over alphabet A . As such, for definite languages, it can be decided whether a given word w belongs to L by looking at the prefix of w of a given fixed length. The same statement holds for *reverse definite* languages, in which we are looking at suffixes. Containing all the definite, or reverse definite, languages is a very weak condition. Note that if a logic can test what is the i -th symbol of a word and is closed under boolean combinations, it already can define all the definite languages. In his paper, Hunt makes an explicit link between intersection emptiness and separability. Hunt writes: “We show that separability is undecidable in general for the same reason that the emptiness-of-intersection problem is undecidable. Therefore, it is unlikely that separability can be used to circumvent the undecidability of the emptiness-of-intersection problem.”

Our Contribution. In this paper, we show that the above mentioned quote does not apply for separability by piecewise testable languages (PTLs): we characterize those full trios for which separability by PTL is decidable. A *full trio* is a nonempty language class that is closed under rational transductions, or, equivalently by Nivat’s theorem (see Nivat, 1968; Berstel, 1979), closed under direct and inverse images under homomorphisms of free monoids, and under intersections with regular languages.

Our characterization states that separability by PTL is decidable if and only if (one of two variants of) an unboundedness problem is decidable. This yields decidability for a range of language classes, such as those with effectively semilinear Parikh images and the languages of labeled vector addition systems. In particular, this means that separability by PTL is decidable for context-free languages, and, according to a very recent result (Hague et al., 2016), even for the languages of higher-order pushdown automata.

The two variants of the unboundedness problem are called *diagonal problem* and *simultaneous unboundedness problem (SUP)*, of which the latter is ostensibly easier than the former. Our reduction of separability by PTL consists of three steps:

- In the first step, we show that (arbitrary) languages I and E are *not* separable by PTL if and only if they possess a certain *common pattern*.
- In the second step we use this fact to reduce the PTL-separation problem to the diagonal problem.
- In the last step, we employ ideal decompositions for downward closed languages to reduce the diagonal problem to the SUP.

For the converse direction, we directly reduce separability by PTL to the SUP.

Another consequence of our characterization is a connection to the problem of *computing downward closures*. It is well known that for any language L , its *downward closure*, i.e., the set of scattered substrings of members of L , is a regular language (Haines, 1969). This is a straightforward consequence of Higman’s Lemma (Higman, 1952). However, given a language L , it is not always possible to compute a finite automaton for the downward closure of L . Since the downward closure appears to be a useful abstraction, this raises the question of when we can compute it. Our characterization here implies that for each full trio \mathcal{S} , downward closures of languages of \mathcal{S} are computable if and only if separability of languages of \mathcal{S} by PTL is decidable.

A curiosity of this work is perhaps the absence of algebraic methods. Most decidability results we are aware of have considered syntactic monoids of regular languages and investigated properties thereof. The exceptions are the recent studies of separability of regular languages by piecewise testable languages (e.g., (Place et al., 2013b; Czerwiński et al., 2013b)).

Here, the situation is different in the sense that for context-free languages the syntactic monoid is infinite and it is difficult to design any algebraic framework for them. So the work shows that it is not always necessary to use algebraic techniques to prove separability questions.

2. PRELIMINARIES AND MAIN RESULTS

The set of all integers and nonnegative integers are denoted by \mathbb{Z} and \mathbb{N} respectively. An *alphabet* is a finite nonempty set, which we usually denote with A , B , or indexed versions thereof. We refer to elements of A as *symbols*. A *word* is a (possibly empty) concatenation $w = a_1 \cdots a_n$ of symbols a_i that come from an alphabet A . A *substring* of w is a sequence $a_j a_{j+1} \cdots a_{j+k}$ of consecutive symbols of w .

The *length* of a word w is n , that is, the number of its symbols. The empty word is denoted by ε . We assume familiarity with regular expressions and regular languages.

The *alphabet of a word* w is the set $\{a_1, \dots, a_n\}$ and is denoted $\text{Alph}(w)$. For a subalphabet $B \subseteq A$, a word $v \in A^*$ is a *B-scattered substring* of w , denoted $v \preceq_B w$, if v can be obtained from w by removing symbols from B , that is, if $v = b_1 \cdots b_m$ and $w \in B^* b_1 B^* \cdots B^* b_m B^*$. Notice that we do not require that $\{b_1, \dots, b_m\} \subseteq B$ or $B \subseteq \{b_1, \dots, b_m\}$. We simply refer to A -scattered substrings as *scattered substrings* and refer to the relation \preceq_A as the *scattered substring relation*, denoted by \preceq . A regular word language over alphabet A is a *piece language* if it is of the form $A^* a_1 A^* \cdots A^* a_n A^*$ for some $a_1, \dots, a_n \in A$, that is, it is the set of words having $a_1 \cdots a_n$ as a scattered substring. A regular language is a *piecewise testable language* if it is a finite Boolean combination of piece languages. The class of all piecewise testable languages is denoted PTL.

2.1. Separability and Common Patterns. The first main result of the paper proves that two (not necessarily regular) languages are not separable by PTL if and only if they have a common pattern. We now make this more precise.

A *factorization pattern* is an element of $(A^*)^{p+1} \times (2^A \setminus \emptyset)^p$ for some $p \geq 0$. In other terms, if (\vec{u}, \vec{B}) is such a factorization pattern, there exist words $u_0, \dots, u_p \in A^*$ and nonempty alphabets $B_1, \dots, B_p \subseteq A$ such that $\vec{u} = (u_0, \dots, u_p)$ and $\vec{B} = (B_1, \dots, B_p)$. For $B \subseteq A$, we denote by B^\circledast the set of words with alphabet exactly B , that is, $B^\circledast = \{w \in B^* \mid \text{Alph}(w) = B\}$. Given a factorization pattern (\vec{u}, \vec{B}) with $\vec{u} = (u_0, \dots, u_p)$ and $\vec{B} = (B_1, \dots, B_p)$, define

$$\mathcal{L}(\vec{u}, \vec{B}, n) = u_0 (B_1^\circledast)^n u_1 \cdots u_{p-1} (B_p^\circledast)^n u_p.$$

In other terms, in a word of $\mathcal{L}(\vec{u}, \vec{B}, n)$, the infix between u_{k-1} and u_k is required to be the concatenation of n words over B_k , each containing *all* symbols of B_k (for each $1 \leq k \leq p$). An infinite sequence $(w_n)_n$ is said to be (\vec{u}, \vec{B}) -adequate if

$$\forall n \in \mathbb{N}, w_n \in \mathcal{L}(\vec{u}, \vec{B}, n).$$

Finally, we say that language L *contains the pattern* (\vec{u}, \vec{B}) if there exists an infinite sequence of words $(w_n)_n$ in L that is (\vec{u}, \vec{B}) -adequate. We can now formally state our first main Theorem which we will prove in Section 3:

Theorem 2.1. *Two word languages I and E are not separable by PTL if and only if they contain a common pattern (\vec{u}, \vec{B}) .*

2.2. Characterizations for Decidable Separability. The second main result is a set of characterizations that say, for *full trios*, when separability by piecewise testable languages is decidable. Full trios, also called *cones*, are language classes that are closed under three operations that we recall next (Berstel, 1979; Ginsburg and Greibach, 1967).

Fix a language L over alphabet A . For an alphabet B , the *B-projection* of a word is its longest scattered substring consisting of symbols from B . The *B-projection* of a language L is the set of all B -projections of words belonging to L . Therefore, the B -projection of L is a language over alphabet $A \cap B$. The *B-upward closure* of a language L is the set of all words that have a B -scattered substring in L , i.e.,

$$\{w \in (A \cup B)^* \mid \exists v \in L \text{ such that } v \preceq_B w\}.$$

In other words, the B -upward closure of L consists of all words that can be obtained by taking a word in L and padding it with symbols from B .

A class of languages \mathcal{C} is *closed* under an operation OP if $L \in \mathcal{C}$ implies that $\text{OP}(L) \in \mathcal{C}$. We use the term *effectively closed* if, furthermore, the representation of $\text{OP}(L)$ can be effectively computed from the representation of L .

A class \mathcal{C} of languages is a *full trio* if it is nonempty and effectively closed under:

- (1) B -projection for every finite alphabet B ,
- (2) B -upward closure for every finite alphabet B , and
- (3) intersection with regular languages.

We note that full trios are usually defined differently (through closures under homomorphisms or rational transductions (Ginsburg and Greibach, 1967; Berstel, 1979)) but we use the above mentioned properties in the proofs and they are easily seen to be equivalent.

We will now introduce three problems whose decidability or computability will be equivalent to decidability of separability by piecewise testable languages: the *diagonal problem*, the *simultaneous unboundedness problem*, and the *downward closure problem*.

Downward Closure Problem. For a language $L \subseteq A^*$, its *downward closure* $L\downarrow$ is defined as the set of all scattered substrings of words in L , that is,

$$L\downarrow = \{u \in A^* \mid \exists v \in L: u \preceq v\}.$$

It is well-known that the downward closure of any language L is regular (Haines, 1969). This is a direct consequence of Higman's Lemma (Higman, 1952), which states that \preceq is a well quasi ordering on A^* , that is, that any infinite sequence of words admits an infinite \preceq -increasing subsequence. Indeed, the complement of a downward closure is closed under A -upward closure. It follows from Higman's Lemma that it has a finite number of \preceq -minimal elements, and that it is actually a finite union of piece languages.

Downward closed languages play an important role in the theory of lossy channel systems (Finkel, 1987; Abdulla and Jonsson, 1996; Finkel and Schnoebelen, 2001), which feature a communication channel where messages can be dropped arbitrarily. These have been investigated extensively, because the WQO property of the subword relation yields decidability results that fail in the non-lossy case (however, the complexity of such decidable problems may be huge (Schnoebelen, 2002; Schmitz and Schnoebelen, 2013)).

Moreover, the downward closure is a useful abstraction of languages: Suppose a language L describes the set of action sequences of a system modeled, for example, by a vector addition system or a pushdown automaton. If this system is observed via a lossy channel, then $L\downarrow$ is the set of sequences seen by the observer (Habermehl et al., 2010). Furthermore, computing a regular representation of $L\downarrow$ for a given language L is in many situations sufficient for safety verification of parametrized asynchronous shared-memory systems (La Torre et al., 2015).

Starting from a regular language given, e.g., by a finite automaton, one can easily compute a representation of its downward closure (one can even obtain precise upper and lower bounds in terms of the size of an automaton recognizing it (Karandikar et al., 2015)). However, despite the fact that the downward closure of a language is always a simple regular language (the complement of a union of piece languages), it is not always possible to effectively compute a finite automaton

for $L\downarrow$ given a description of L . Such negative results are known, for example, for Church-Rosser languages (Gruber et al., 2007) and reachability sets of lossy channel systems (Mayr, 2003). We say that *downward closures are computable for \mathcal{C}* if a finite automaton for $L\downarrow$ can be algorithmically computed for every language L in \mathcal{C} .

Unboundedness Problems. Let $A = \{a_1, \dots, a_n\}$ ordered with $a_1 < \dots < a_n$. For a symbol $a \in A$ and a word $w \in A^*$, let $\#_a(w)$ denote the number of occurrences of a in w . The *Parikh image* of a word w is the n -tuple $(\#_{a_1}(w), \dots, \#_{a_n}(w))$. The *Parikh image* of a language L is the set of all Parikh images of words from L . A tuple $(m_1, \dots, m_n) \in \mathbb{N}^n$ is *dominated* by a tuple $(d_1, \dots, d_n) \in \mathbb{N}^n$ if $d_i \geq m_i$ for every $i = 1, \dots, n$.

Definition 2.2. *The diagonal problem for \mathcal{C} is the following decision problem:*

Input: A language $L \subseteq A^*$ from \mathcal{C} .

Question: Is each tuple $(m, \dots, m) \in \mathbb{N}^n$ dominated by some tuple in the Parikh image of L ?

Equivalently, the diagonal problem for \mathcal{C} asks whether there are *infinitely many* tuples (m, \dots, m) dominated by some tuple in the Parikh image of L .

The *simultaneous unboundedness problem (SUP)* is a restricted version of the diagonal problem where the input is a language in which the symbols are grouped together:

Definition 2.3. *The simultaneous unboundedness problem (SUP) for \mathcal{C} is the following decision problem:*

Input: A language $L \subseteq b_1^* \dots b_n^*$ from \mathcal{C} , for some ordering b_1, \dots, b_n of A .

Question: Is each tuple $(m, \dots, m) \in \mathbb{N}^n$ dominated by some tuple in the Parikh image of L ?

In other words, the SUP asks whether $L\downarrow = b_1^* \dots b_n^*$.

Characterizations. Perhaps surprisingly, our second main result here implies that in a full trio, separability by PTL is decidable if and only if downward closures are computable. Our proof relies on the following characterization of computability of downward closures:

Theorem 2.4 (Zetsche, 2015a). *Let \mathcal{C} be a full trio. Then downward closures are computable for \mathcal{C} if and only if the SUP is decidable for \mathcal{C} .*

We are now ready to state the second main result, which is an extension of Theorem 2.4 and connects it to separability:

Theorem 2.5. *For each full trio \mathcal{C} , the following are equivalent:*

- (1) *Separability of \mathcal{C} by PTL is decidable.*
- (2) *The diagonal problem for \mathcal{C} is decidable.*
- (3) *The SUP for \mathcal{C} is decidable.*
- (4) *Downward closures are computable for \mathcal{C} .*

Theorem 2.4 provides the equivalence between 3 and 4. We prove the remaining equivalences in Section 4. In particular, we present an algorithm to decide separability for full trios that have a decidable diagonal problem, showing one direction of the equivalence. The algorithm does not rely on semilinearity of Parikh images. For example, in Section 5 we apply the theorem to Vector Addition System languages, which do not have a semilinear Parikh image.

3. COMMON PATTERNS

In this section we prove Theorem 2.1. We say that an infinite sequence is *adequate* if it is (\vec{u}, \vec{B}) -adequate for some factorization pattern. The following statement can be shown using Simon's Factorization Forest Theorem (Simon, 1990).

Lemma 3.1. *Every infinite sequence $(w_n)_n$ of words admits an adequate subsequence.*

Proof. We use Simon's Factorization Forest Theorem, which we recall. See (Simon, 1990; Kufleitner, 2008; Bojańczyk, 2009; Colcombet, 2010, 2016) for proofs and extensions of this theorem. A *factorization tree* of a nonempty word x is a finite ordered unranked tree $T(x)$ whose nodes are labeled by nonempty words, and such that:

- all leaves of $T(x)$ are labeled by symbols,
- all internal nodes of $T(x)$ have at least 2 children, and
- if a node labeled y has k children labeled y_1, \dots, y_k from left to right, then $y = y_1 \cdots y_k$.

Given a semigroup morphism $\varphi : A^+ \rightarrow S$ into a finite semigroup S , such a factorization tree is called φ -*Ramseyan* if every internal node has either 2 children, or k children labeled y_1, \dots, y_k , in which case φ maps all words y_1, \dots, y_k to the same idempotent of S , i.e., to an element e such that $ee = e$. Simon's Factorization Forest Theorem states that every word has a φ -Ramseyan factorization tree of height at most $3|S|$.

Let $(w_n)_n$ be an infinite sequence of words. We use Simon's Factorization Forest Theorem with the morphism $\text{Alph} : A^+ \rightarrow 2^A$. Recall that Alph maps a word w to the set of symbols used in w .

Consider a sequence $(T(w_n))_n$, where $T(w_n)$ is an Alph -Ramseyan tree of w_n , given by the Factorization Forest Theorem. In particular, one may choose $T(w_n)$ of depth at most $3 \cdot 2^{|A|}$. Therefore, extracting a subsequence if necessary, one may assume that the sequence of depths of the trees $T(w_n)$ is a constant H . We argue by induction on H . If $H = 0$, then every w_n is a symbol. Hence, one may extract from $(w_n)_n$ a constant subsequence, say $(w)_{n \in \mathbb{N}}$, which is therefore $((w), ())$ -adequate, that is, adequate for the factorization pattern consisting of (w) and the empty vector $()$. Hence, it is adequate.

We denote the arity of the root of $T(w_n)$ by $\text{arity}(w_n)$ and we call it the arity of w_n . If $H > 0$, two cases may arise:

- (1) One can extract from $(w_n)_n$ a subsequence of bounded arity. Therefore, one may extract a subsequence of constant arity, say K , from w_n . This implies that each w_n can be written as a concatenation of K words

$$w_n = w_{n,1} \cdots w_{n,K},$$

where $w_{n,i}$ is the label of the i -th child of the root in $T(w_n)$. Therefore, the Alph -Ramseyan subtree of each $w_{n,i}$ is of height at most $H - 1$. By induction, one can extract from every $(w_{n,i})_n$ an adequate subsequence. Proceeding iteratively for $i = 1, 2, \dots, K$, one extracts from $(w_n)_n$ a subsequence $(w_{\sigma(n)})_n$ such that every $(w_{\sigma(n),i})_n$ is adequate. But a finite concatenation of adequate sequences is obviously adequate. Therefore, the subsequence $(w_{\sigma(n)})_n$ of $(w_n)_n$ is also adequate.

- (2) The arity of w_n grows to infinity. Therefore, extracting if necessary, one can assume for every n , that $\text{arity}(w_n) \geq \max(n, 3)$. Since each factorization tree is Alph -Ramseyan and since the arity of the root of each tree is at least 3, all children of the root of the n -th tree map to the same idempotent in 2^A , say $B_n \subseteq A$. Since 2^A is finite, one can further extract a subsequence, say $w_{\sigma(n)}$, such that $B_{\sigma(n)}$ is constant, equal to some $B \subseteq A$. To sum up, each word from the subsequence is of the form

$$w_{\sigma(n)} = w_{n,1} \cdots w_{n,K_n},$$

with $K_n \geq n$ and, where the alphabet of $w_{n,i}$ is B . Therefore, $w_{\sigma(n)} \in (B^*)^{K_n} \subseteq (B^*)^n$. Therefore, $(w_{\sigma(n)})_n$ is $((\varepsilon, \varepsilon), (B))$ -adequate, hence it is adequate. \square

For a word w , denote its first (resp., last) symbol by $\text{first}(w)$, (resp., $\text{last}(w)$). We call a factorization pattern $(\vec{u}, \vec{B}) = ((u_0, \dots, u_p), (B_1, \dots, B_p))$ *proper* if

- (1) for all $i = 0, \dots, p-1$, we have $\text{last}(u_i) \notin B_{i+1}$, for all $i = 1, \dots, p$, we have $\text{first}(u_i) \notin B_i$, and
- (2) for all $i = 1, \dots, p-1$, if $u_i = \varepsilon$, then we have $(B_i \not\subseteq B_{i+1} \text{ and } B_{i+1} \not\subseteq B_i)$.

Note that if a sequence $(w_n)_n$ is adequate, then there exists a proper factorization pattern (\vec{u}, \vec{B}) such that $(w_n)_n$ is (\vec{u}, \vec{B}) -adequate. This is easily seen from the following observations and their symmetric counterparts:

$$\begin{aligned} u = a_1 \cdots a_k \text{ and } a_k \in B &\Rightarrow a_1 \cdots a_k (B^\otimes)^n \subseteq a_1 \cdots a_{k-1} (B^\otimes)^n, \\ B_i \subseteq B_{i+1} &\Rightarrow (B_i^\otimes)^n (B_{i+1}^\otimes)^n \subseteq (B_{i+1}^\otimes)^n. \end{aligned}$$

The following lemma gives a condition under which two sequences share a factorization pattern and is very similar to (Almeida, 1994, Theorem 8.2.6). In its statement, we write $v \sim_n w$ for two words v and w if they have the same scattered substrings up to length n , that is, for every word u of length at most n , we have $u \preceq v$ if and only if $u \preceq w$. Notice that \sim_n is an equivalence relation for every $n \in \mathbb{N}$.

Lemma 3.2. *Let (\vec{u}, \vec{B}) and (\vec{t}, \vec{C}) be proper factorization patterns. Let $(v_n)_n$ and $(w_n)_n$ be two sequences of words such that*

- $(v_n)_n$ is (\vec{u}, \vec{B}) -adequate
- $(w_n)_n$ is (\vec{t}, \vec{C}) -adequate, and
- $v_n \sim_n w_n$ for every $n \geq 0$.

Then, $\vec{u} = \vec{t}$ and $\vec{B} = \vec{C}$.

Proof. For a factorization pattern (\vec{u}, \vec{B}) , we define

$$\|(\vec{u}, \vec{B})\| = \left(\sum_{i=0}^{\ell} |u_i| \right) + \ell,$$

where $\ell + 1$ is the length of the vector $\vec{u} = (u_0, \dots, u_\ell)$. Let

$$k = \max(\|(\vec{u}, \vec{B})\|, \|(\vec{t}, \vec{C})\|) + 1.$$

Consider the first word of the sequence $(v_n)_n$, i.e., $v_0 = u_0 r_1 u_1 \cdots r_\ell u_\ell$, where $\text{Alph}(r_i) = B_i$. Define

$$(1) \quad v_0^{(k)} = u_0 r_1^k u_1 \cdots r_\ell^k u_\ell.$$

Recall that $(v_n)_n$ being a (\vec{u}, \vec{B}) -adequate sequence means that

$$(2) \quad \forall n \in \mathbb{N}, \quad v_n \in u_0 (B_1^\otimes)^n u_1 \cdots u_{\ell-1} (B_\ell^\otimes)^n u_\ell.$$

Let $N = k \cdot \max(|r_1|, \dots, |r_n|)$. By (1) and (2), we get $v_0^{(k)} \preceq v_N$. Let $M \geq \max(N, |v_0^{(k)}|)$, so that $v_0^{(k)}$ is a scattered substring of v_M of length at most M . Since $v_M \sim_M w_M$, this gives that

$$(3) \quad v_0^{(k)} \preceq w_M.$$

To show that $(\vec{u}, \vec{B}) = (\vec{t}, \vec{C})$, we first define a bijection between elements of the sequence of indexed alphabets in \vec{B} and elements of the sequence of those in \vec{C} . For this, we embed $v_0^{(k)}$ in w_M in the ‘leftmost’ way. Let us make this precise: let $x, y \in A^*$ with $x = x_1 \cdots x_{|x|}$ and $y = y_1 \cdots y_{|y|}$, where $x_i, y_i \in A$. If $x \preceq y$, by definition there exists an increasing mapping $h : \{1, \dots, |x|\} \rightarrow \{1, \dots, |y|\}$ such that $x_i = y_{h(i)}$ holds for all i . We call such a mapping an *embedding* of x in y . The *leftmost embedding* h_{left} of x in y is such that additionally, for any embedding h , we have $h_{\text{left}}(i) \leq h(i)$ for all i .

To define the bijection, set $\mathbf{B} = \{(B_1, 1), \dots, (B_\ell, \ell)\}$ and $\mathbf{C} = \{(C_1, 1), \dots, (C_m, m)\}$, where m denotes the length of the vector \vec{C} . We define a function $f : \mathbf{B} \rightarrow \mathbf{C}$ as follows. Consider the leftmost embedding h_{left} of $v_0^{(k)}$ in w_M , which exists by (3). Since $(w_n)_n$ is a (\vec{t}, \vec{C}) -adequate sequence, one may write

$$w_M = t_0 s_1 t_1 \cdots t_{m-1} s_m t_m \quad \text{with } s_j \in (C_j^*)^M \text{ for all } j.$$

For i in $\{1, \dots, \ell\}$, consider the substring r_i^k of $v_0^{(k)}$. By definition of $k > \|(\vec{t}, \vec{C})\|$ and since $|r_i| > 0$, the pigeonhole principle implies that all positions of one of these k copies of r_i must be mapped by h_{left} to positions of some substring $s_j \in (C_j^*)^M$. We define $f(B_i, i) = (C_j, j)$ for the smallest such index j .

The function $g : \mathbf{C} \rightarrow \mathbf{B}$ is defined symmetrically, by exchanging the roles of $(v_n)_n$ and $(w_n)_n$. Observe that the functions f and g preserve the order of the indices. Moreover, since $\text{Alph}(r_i) = B_i$ and $r_i \preceq s_j \in (C_j^*)^M$, we obtain that $f(B_i, i) = (C_j, j)$ entails $B_i \subseteq C_j$. Similarly, if $g(C_j, j) = (B_i, i)$, then $C_j \subseteq B_i$. If we show that f and g define a bijective correspondence between \mathbf{B} and \mathbf{C} , then $\ell = m$. The above observations would then imply that $B_i = C_i$, for every i .

To establish that f and g are inverses of one another, we apply Lemma 8.2.5 from (Almeida, 1994), which we shall first repeat:

Lemma 3.3 (Almeida, 1994, Lemma 8.2.5). *Let X and Y be finite sets and let P be a partially ordered set. Let $f : X \rightarrow Y, g : Y \rightarrow X, p : X \rightarrow P$ and $q : Y \rightarrow P$ be functions such that*

- (1) *for any $x \in X, p(x) \leq q(f(x))$,*
- (2) *for any $y \in Y, q(y) \leq p(g(y))$,*
- (3) *if $x_1, x_2 \in X, f(x_1) = f(x_2)$ and $p(x_1) = q(f(x_1))$, then $x_1 = x_2$,*
- (4) *if $y_1, y_2 \in Y, g(y_1) = g(y_2)$ and $q(y_1) = p(g(y_1))$, then $y_1 = y_2$.*

Then f and g are mutually inverse functions and $p = q \circ f$ and $q = p \circ g$.

To apply this lemma, set $X = \mathbf{B}, Y = \mathbf{C}, P$ as the set of subalphabets of A partially ordered by inclusion, and let p and q be the projections onto the first coordinate.

Let us verify that f and g fulfill all conditions of Lemma 3.3. Item 1 holds since for all i and j , $f(B_i, i) = (C_j, j)$ implies that $B_i \subseteq C_j$. Item 2 holds symmetrically.

For Item 3, suppose that $f(B_{i_1}, i_1) = f(B_{i_2}, i_2) = (C_j, j)$ and that $p(B_{i_1}, i_1) = q(f(B_{i_1}, i_1))$, that is, $B_{i_1} = C_j$. The first condition, $f(B_{i_1}, i_1) = f(B_{i_2}, i_2) = (C_j, j)$, means that a substring r_{i_1} and a substring r_{i_2} of $v_0^{(k)}$ have all their positions mapped by h_{left} to positions of s_j in w_M . Therefore, all intermediate positions are also mapped to positions of s_j , which implies, using the second condition, $B_{i_1} = C_j$, that $\text{Alph}(r_{i_1} u_{i_1} \cdots r_{i_2}) \subseteq \text{Alph}(s_j) = C_j = B_{i_1} = \text{Alph}(r_{i_1})$. But we assumed that (\vec{u}, \vec{B}) is a *proper* factorization pattern, so i_1 must be equal to i_2 . This shows that item 3 holds. Item 4 is dual.

It follows that indeed f and g define a bijective correspondence between \mathbf{B} and \mathbf{C} , thus $\ell = m$ and $B_i = C_i$, for every i . Since we are dealing with proper factorization patterns, $v_0^{(k)} \preceq w_M$ now implies that $t_i \preceq u_i$, for every i . Symmetrically, $u_i \preceq t_i$ for every i . Thus, for every i , $u_i = t_i$. \square

Proof of Theorem 2.1. *Two word languages I and E are not separable by PTL if and only if they contain a common pattern (\vec{u}, \vec{B}) .*

Proof. We rely on the following characterization of PTL-separation: I and E are *not* PTL-separable if and only if for every $n \in \mathbb{N}$, there exist $v_n \in I$ and $w_n \in E$ such that $v_n \sim_n w_n$. Indeed, suppose that there exists $n \in \mathbb{N}$ such that there are no words in I (respectively, E) that are \sim_n -equivalent. Then $[I]_n := \{v \mid \exists w \in I \text{ such that } v \sim_n w\}$ separates I from E . It remains to verify that it is a PTL language. Observe that \sim_n has finite index, because a \sim_n -class is defined by a subset of words

of length at most n . Therefore, $[I]_n$ is a finite union of \sim_n -classes. To show that $[I]_n$ is PTL, it is therefore enough to check that a single \sim_n -class is PTL, which is the case since the \sim_n -class of u is

$$\bigcap_{v \preceq u, |v| \leq n} v\uparrow \cap \bigcap_{v \not\preceq u, |v| \leq n} A^* \setminus (v\uparrow)$$

where $v\uparrow$ is the A -upward closure of $\{v\}$, i.e., the piece language $A^*a_1A^* \cdots A^*a_kA^*$, if $v = a_1 \cdots a_k$.

On the other hand, suppose that for every $n \in \mathbb{N}$, there are words in I (respectively, E) that are \sim_n -equivalent. A piecewise testable language is a union of \sim_n -equivalence classes for some n , therefore any piecewise testable language containing I will, in this case, have non-empty intersection with E .

To prove the “if”-direction, we show that for every n , there exist words in I and in E that are \sim_n -equivalent. By hypothesis, I and E contain a common factorization pattern (\vec{u}, \vec{B}) . By definition, this gives us two sequences of words $(v_n)_n$ and $(w_n)_n$ such that for all n ,

$$\begin{aligned} v_n &\in I \cap u_0(B_1^\otimes)^n u_1 \cdots u_{p-1}(B_p^\otimes)^n u_p, \\ w_n &\in E \cap u_0(B_1^\otimes)^n u_1 \cdots u_{p-1}(B_p^\otimes)^n u_p. \end{aligned}$$

Observe that for all i , $(B_i^\otimes)^n$ contains precisely all words from $B_i^{\leq n}$ as scattered substrings of size up to n . It follows that $v_n \sim_n w_n$.

We now prove the “only-if” direction. The existence of $v_n \sim_n w_n$ for every $n \in \mathbb{N}$ defines an infinite sequence of pairs $(v_n, w_n)_n$, from which we will iteratively extract infinite subsequences to obtain additional properties, while keeping \sim_n -equivalence. By Lemma 3.1, one can extract from $(v_n, w_n)_n$ a subsequence whose first component forms an adequate sequence. From this subsequence of pairs, using Lemma 3.1 again, we extract a subsequence whose second component is also adequate (note that the first component remains adequate). Therefore, one can assume that both $(v_n)_n$ and $(w_n)_n$ are themselves adequate. This means there exist proper factorization patterns for which $(v_n)_n$ resp. $(w_n)_n$ are adequate. Lemma 3.2 shows that one can choose the *same* proper factorization pattern (\vec{u}, \vec{B}) such that both $(v_n)_n$ and $(w_n)_n$ are (\vec{u}, \vec{B}) -adequate. This means that I and E contain a common pattern (\vec{u}, \vec{B}) . \square

4. THE CHARACTERIZATION FOR SEPARABILITY

In this section we prove our main characterization:

Theorem 2.5. *For each full trio \mathcal{C} , the following are equivalent:*

- (1) *Separability of \mathcal{C} by PTL is decidable.*
- (2) *The diagonal problem for \mathcal{C} is decidable.*
- (3) *The SUP for \mathcal{C} is decidable.*
- (4) *Downward closures are computable for \mathcal{C} .*

The equivalence between 3 and 4 is immediate from Theorem 2.4. The implication “2 \Rightarrow 3” is trivial because the SUP is a special case of the diagonal problem. We prove the other direction “3 \Rightarrow 2” and the implication “1 \Rightarrow 3” in Section 4.1. Finally, we prove the implication “2 \Rightarrow 1” in Section 4.2, by giving an algorithm for separability if the diagonal problem is decidable.

4.1. Algorithms for the Diagonal Problem and the SUP. In this section, we reduce the diagonal problem to the simultaneous unboundedness problem and to separability by PTL. It should be noted that the implication “3 \Rightarrow 2” already follows easily from the equivalence between 3 and 4 (shown in (Zetzsche, 2015a)): when one can compute downward closures, the diagonal problem reduces to the case of regular languages. However, the algorithm for downward closure computation in (Zetzsche, 2015a) is not accompanied by any complexity bounds. Therefore, we mention here a reduction that can be carried out in non-deterministic polynomial time if we assume that the full

trio operations require only polynomial time (Lemma 4.2). This reduction is based on the fact that downward closed languages can be written as finite unions of ideals, which is also a central ingredient in (Zetsche, 2015a). An *ideal* is a language of the form $B_0^* \{b_1, \varepsilon\} B_1^* \cdots \{b_m, \varepsilon\} B_m^*$, where $m \geq 0$, b_1, \dots, b_m are symbols, and B_0, \dots, B_m are alphabets. Clearly, every ideal is downward closed. We use the following result by Jullien, which has later been rediscovered in (Abdulla et al., 2004).

Theorem 4.1 (Jullien, 1969). *Every downward closed language is a finite union of ideals.*

In fact, there is a notion of ideal in well-quasi-orderings for which this result holds in general. See for instance (Finkel and Goubault-Larrecq, 2009; Leroux and Schmitz, 2015) for other perspectives. We prove the implication “3 \Rightarrow 2” of Theorem 2.5.

Lemma 4.2. *Let \mathcal{C} be a full trio. If the SUP is decidable for \mathcal{C} , then the diagonal problem is decidable for \mathcal{C} as well.*

Proof. Let $L \subseteq A^*$ with $A = \{a_1, \dots, a_n\}$. We say that L satisfies the *diagonal property* if for each $m \in \mathbb{N}$, the vector (m, \dots, m) (with $|A|$ entries) is dominated by some vector in the Parikh image of L . We claim that L satisfies the diagonal property if and only if we can order the symbols of A as $A = \{b_1 \leq \dots \leq b_n\}$ such that $b_1^* \cdots b_n^* \subseteq L\downarrow$. The “if” direction is immediate, so suppose L satisfies the diagonal property.

Since L satisfies the diagonal property, so does $L\downarrow$. According to Theorem 4.1, we can write $L\downarrow$ as a finite union of ideals. This means that at least one of these ideals has to satisfy the diagonal property. Let $I = C_0^* \{c_1, \varepsilon\} C_1^* \cdots \{c_m, \varepsilon\} C_m^*$ be one such ideal. Then we have $C_0 \cup \dots \cup C_m = A$, since any element of $A \setminus (C_0 \cup \dots \cup C_m)$ can occur at most m times in words in I . Hence, if we order the elements of A linearly by picking first the elements of C_0 , then those of $C_1 \setminus C_0$, and so forth, the resulting ordering $A = \{b_1 \leq \dots \leq b_n\}$ clearly satisfies $b_1^* \cdots b_n^* \subseteq I \subseteq L\downarrow$. This proves our claim.

Assume that \mathcal{C} is a full trio for which the SUP is decidable. We show that the diagonal problem is decidable for \mathcal{C} as well. Let $L \subseteq A^*$ be a language. We guess a linear ordering $A = \{b_1 \leq \dots \leq b_n\}$ and construct the language $K = L\downarrow \cap b_1^* \cdots b_n^*$. Notice that $L\downarrow$ and therefore K can be constructed due to Theorem 2.4. Then, K is an instance of the SUP. By our claim, L satisfies the diagonal property if and only if there exists a linear ordering $A = \{b_1 \leq \dots \leq b_n\}$ such that $K = L\downarrow \cap b_1^* \cdots b_n^*$ is a positive instance of the SUP. \square

This concludes the proof of “3 \Rightarrow 2”, of Theorem 2.5, so that we now have established that Items 2, 3 and 4 are equivalent. It remains to connect these problems with the separation problem.

The correctness proof of our reduction of the SUP to the separability problem employs a characterization of separability by PTL from (Czerwiński et al., 2013b). For languages $K, L \subseteq A^*$, a (K, L) -zigzag is an infinite sequence of words $(w_i)_{i \in \mathbb{N}}$ such that

- (i) $w_i \preceq w_{i+1}$ for every $i \in \mathbb{N}$,
- (ii) $w_i \in K$ for every even i and
- (iii) $w_i \in L$ for every odd i .

The characterization for separability by PTL is then as follows.

Theorem 4.3 (Czerwiński et al., 2013b). *Let $K, L \subseteq A^*$ be languages. Then L and K are separable by PTL if and only if there is no (K, L) -zigzag.*

In the following lemma, we use the definition of full trios as being closed under rational transductions. A subset $T \subseteq A^* \times B^*$ for alphabets A, B is called *rational transduction* if there is an alphabet C , a regular language $R \subseteq C^*$, and homomorphisms $\alpha: C^* \rightarrow A^*$, $\beta: C^* \rightarrow B^*$ such that we have $T = \{(\alpha(w), \beta(w)) \mid w \in R\}$. For a language $L \subseteq A^*$ and a rational transduction

$T \subseteq A^* \times B^*$, we define

$$TL = \{w \in B^* \mid \exists v \in L: (v, w) \in T\}.$$

It is well-known that a language class \mathcal{C} is a full trio if and only if for each $L \in \mathcal{C}$ and each rational transduction T , we have effectively $TL \in \mathcal{C}$, see (Berstel, 1979) for instance.

Lemma 4.4. *Let \mathcal{C} be a full trio. If separability by PTL is decidable for \mathcal{C} , then the SUP is decidable for \mathcal{C} as well.*

Proof. Let $A = \{a_1, \dots, a_n\}$ and let $L \subseteq a_1^* \cdots a_n^*$ be a language in \mathcal{C} . Let $T \subseteq A^* \times A^*$ be the rational transduction

$$T = \{(a_1^{k_1} \cdots a_n^{k_n}, a_1^{2k_1} \cdots a_n^{2k_n}) \mid k_1, \dots, k_n \geq 0\}.$$

Moreover, consider the regular language

$$K = \{a_1^{2k_1+1} \cdots a_n^{2k_n+1} \mid k_1, \dots, k_n \geq 0\}.$$

We claim that $T(L\downarrow)$ and K are inseparable by PTL if and only if $L\downarrow = a_1^* \cdots a_n^*$. We first show that this claim implies the lemma. Indeed, since \mathcal{C} is a full trio, we can effectively compute representations (in the form of a finite automaton or in the form customary for \mathcal{C}) of $T(L\downarrow)$. Since K is a member of \mathcal{C} (since every full trio contains the family of regular languages), the claim clearly implies the lemma.

We now prove the claim. Suppose that $L\downarrow = a_1^* \cdots a_n^*$. Then we have $T(L\downarrow) = (a_1 a_1)^* \cdots (a_n a_n)^*$ and the sequence $(w_i)_{i \in \mathbb{N}}$ with $w_i = a_1^i \cdots a_n^i$ is a $(T(L\downarrow), K)$ -zigzag, meaning $T(L\downarrow)$ and K are inseparable by PTL, by Theorem 4.3.

Now suppose $T(L\downarrow)$ and K are inseparable by PTL. Then again by Theorem 4.3, there is a $(T(L\downarrow), K)$ -zigzag $(w_i)_{i \in \mathbb{N}}$. This means, for $i \in \{1, \dots, n\}$, we have $w_i \preceq w_{i+1}$, so that $|w_i|_a \leq |w_{i+1}|_a$ for $a \in A$. By construction of T and K , the numbers $|w_i|_a$ and $|w_{i+1}|_a$ are incongruent modulo 2, which implies $|w_i|_a < |w_{i+1}|_a$ and therefore $|w_i|_a \geq i$. Moreover, since the sequence is a zigzag, we have $\{w_{2i} \mid i \geq 0\} \subseteq T(L\downarrow)$ and thus $L\downarrow = a_1^* \cdots a_n^*$. This completes the proof of our lemma. \square

4.2. The Algorithm for Separability. It only remains to prove the implication “ $2 \Rightarrow 1$ ” to finish the proof of Theorem 2.5. In this section we show that, for full trios with decidable diagonal problem, we can decide separability by PTL. Fix two languages I and E from a full trio \mathcal{C} which has decidable diagonal problem.

To test whether I is separable from E by a piecewise testable language S , we run two semi-procedures in parallel. The *positive* one looks for a witness that I and E are separable by PTL, whereas the *negative* one looks for a witness that they are *not* separable by a PTL. Since one of the semi-procedures always terminates, we have an effective algorithm that decides separability. It remains to describe the two semi-procedures.

Positive semi-procedure. We first note that, when a full trio has decidable diagonal problem, it also has decidable emptiness. Indeed, emptiness of a language $L \subseteq A^*$ can be decided by taking the A -upward closure of L (which can be effectively computed from L , since it can be implemented by a rational transduction). In the resulting language, the diagonal problem returns true if and only if L is nonempty.

The positive semi-procedure enumerates all PTLs over the union of the alphabets of I and E . For every PTL S it checks whether S is a separator, so if $I \subseteq S$ and $E \cap S = \emptyset$. The first test is equivalent to $I \cap (A^* \setminus S) = \emptyset$. Thus both tests boil down to checking whether the intersection of a language from the class \mathcal{C} (I or E , respectively) and a regular language (S and $A^* \setminus S$, respectively) is empty. This is decidable, as \mathcal{C} is effectively closed under taking intersections with regular languages and has decidable emptiness problem.

Negative semi-procedure. Theorem 2.1 shows that there is always a finite witness for inseparability: a pattern (\vec{u}, \vec{B}) . The negative semi-procedure enumerates all possible patterns and for each one, checks the condition of Theorem 2.1. We now show how to test this condition, i.e., for a pattern (\vec{u}, \vec{B}) test whether for all $n \in \mathbb{N}$ the intersection of $L(\vec{u}, \vec{B}, n)$ with both I and E is nonempty.

Checking the condition. Here we show for an arbitrary language from \mathcal{C} how to check whether for all $n \in \mathbb{N}$ its intersection with the language $L(\vec{u}, \vec{B}, n)$ is nonempty. Fix $L \in \mathcal{C}$ over an alphabet A and a pattern (\vec{u}, \vec{B}) , where $\vec{u} = (u_0, \dots, u_k)$ and $\vec{B} = (B_1, \dots, B_k)$. Intuitively, we just consider a diagonal problem with some artifacts: we are counting the number of occurrences of alphabets B_i and checking whether those numbers can simultaneously become arbitrarily big.

We show decidability of the non-separability problem by a formal reduction to the diagonal problem. We perform a sequence of steps. In every step we will slightly modify the considered language L and appropriately customize the condition to be checked. Using the closure properties of the class \mathcal{C} we will ensure that the investigated language still belongs to \mathcal{C} .

First we add special symbols $\$i$, for $i \in \{1, \dots, k\}$, which do not occur in A . These symbols are meant to count how many times alphabet B_i is “fully occurring” in the word. Then we will assure that words are of the form

$$u_0 (B_1 \cup \{\$1\})^* u_1 \cdots u_{k-1} (B_k \cup \{\$k\})^* u_k,$$

which already is close to what we need for the pattern. Then we will check that between every two symbols $\$i$ (with the same i), every symbol from B_i occurs, so that the $\$i$ are indeed counting the number of iterations through the entire alphabet B_i . Finally we will remove all the symbols except those from $\{\$1, \dots, \$k\}$. The resulting language will contain only words of the form $\$1^* \$2^* \cdots \$k^*$ and the condition to be checked will be exactly the diagonal problem.

More formally, let $L_0 := L$. We modify iteratively L_0 , resulting in L_1, L_2, L_3 , and L_4 . Each of them will be in \mathcal{C} and we describe them next.

Language L_1 is the $\{\$1, \dots, \$k\}$ -upward closure of L_0 . Thus, L_1 contains, in particular, all words where the $\$i$ are placed “correctly”, i.e., in between two $\$i$ -symbols the whole alphabet B_i should occur. However at this moment we do not check it. By closure under B -upward closures, which can be implemented by rational transductions, language L_1 belongs to \mathcal{C} .

Note that L_1 also contains words in which the $\$i$ -symbols are placed totally arbitrary. In particular, they can occur in the wrong order. The idea behind L_2 is to consider only those words in which the $\$i$ -symbols were placed at least in the right areas. Concretely, L_2 is the intersection of L_1 with the language

$$u_0 (B_1 \cup \{\$1\})^* u_1 \cdots u_{k-1} (B_k \cup \{\$k\})^* u_k.$$

Since \mathcal{C} is a full trio, it is closed under intersection with regular languages, whence L_2 belongs to \mathcal{C} .

Language L_2 may still contain words, such that in between two $\$i$ -symbols not *all* the symbols from B_i occur. We get rid of these by intersecting L_2 with the regular language

$$u_0 (\$1 B_1^*)^* \$1 u_1 \cdots u_{k-1} (\$k B_k^*)^* \$k u_k.$$

As such, we obtain L_3 which, again by closure under intersection with regular languages, belongs to \mathcal{C} .¹

Note that the intersection of $L = L_0$ with the language $L(\vec{u}, \vec{B}, n)$ is nonempty if and only if L_3 contains a word with precisely $n+1$ symbols $\$i$ for every $i \in \{1, \dots, k\}$. Indeed, L_3 just contains the (slightly modified versions of) words from L_0 which fit into the pattern and in which the symbols $\$i$ “count” occurrences of B_i^* . Furthermore, for every word in L_3 , the word obtained by removing

¹Of course, one could also immediately obtain L_3 from L_1 by performing a single intersection with a regular language.

some occurrences of some $\$i$ is in L_3 as well. It is thus enough to focus on the $\$i$ -symbols. Language L_4 is therefore the $\{\$, \dots, \$k\}$ -projection of L_3 . By the closure under projections, language L_4 belongs to \mathcal{C} . The words contained in L_4 are therefore of the form

$$\$_1^{a_1} \dots \$_k^{a_k},$$

such that there exists $w \in L$ with at least $a_i - 1$ occurrences of substrings of B_i^\otimes between u_{i-1} and u_i . Therefore, $L \cap L(\vec{u}, \vec{B}, n)$ is nonempty for all $n \geq 0$ if and only if the tuple (n, \dots, n) belongs to the Parikh image of L_4 for infinitely many $n \geq 0$. This is precisely the diagonal problem, which we know to be decidable for \mathcal{C} . \square

5. DECIDABLE CLASSES

In this section we show that separability by piecewise testable languages is decidable for a wide range of classes, by proving that they meet the conditions of Theorem 2.5, in particular, for context-free languages and languages of labeled vector addition systems (which are the same as languages of labeled Petri nets).

Effectively semilinear Parikh images. A number of language classes is known to exhibit effectively semilinear Parikh images. A set $S \subseteq \mathbb{N}^k$ is *linear* if it is of the form

$$S = \{v + n_1 v_1 + \dots + n_m v_m \mid n_1, \dots, n_m \in \mathbb{N}\}$$

for some *base* vector $v \in \mathbb{N}^k$ and *period* vectors $v_1, \dots, v_m \in \mathbb{N}^k$. A *semilinear* set is a finite union of linear sets. We say that a full trio \mathcal{C} exhibits *effectively semilinear Parikh images* if every language in \mathcal{C} has a semilinear Parikh image and one can compute a representation as a (finite) union of linear sets.

Clearly, one can decide the diagonal problem for language classes with effectively semilinear Parikh images. This amounts to checking whether in a representation of the Parikh image of the given language, there is some linear set in which each of the symbols occurs in some period vector.

Another option is to use Presburger logic. Semilinear sets are exactly those definable by Presburger logic. Moreover, the translation is effective. Assume that $|A| = k$, so the Parikh image P of the considered language is a subset of \mathbb{N}^k and ϕ is a Presburger formula describing P having k free variables. Then

$$\psi = \forall_{n \in \mathbb{N}} \exists_{x_1, x_2, \dots, x_k} \left(\bigwedge_{i \in \{1, \dots, k\}} (x_i \geq n) \right) \wedge \phi(x_1, x_2, \dots, x_k)$$

is true if and only if the diagonal problem for the considered language is answered positively. Decidability of the Presburger logic finishes the proof of decidability of the diagonal problem. We refer for the details of semilinear sets and Presburger logic to (Ginsburg and Spanier, 1966).

Examples of full trios with effectively semilinear Parikh images are *context-free languages* (Parikh, 1966), *multiple context-free languages* (Seki et al., 1991), languages of *reversal-bounded counter automata* (Ibarra, 1978), *stacked counter automata* (Zetzsche, 2015b), and *finite index matrix languages* Dassow and Păun (1989).

Higher-Order Pushdown Automata. Very recently, it was shown that the diagonal problem is decidable for higher-order pushdown automata (Hague et al., 2016). The authors use an inductive approach: The diagonal problem for order- $(n+1)$ pushdown automata is decided by constructing an order- n pushdown automaton that is equivalent with respect to the diagonal problem. Hence, the algorithm arrives at an order-1 pushdown automata, which is an ordinary pushdown automaton.

Languages of Labeled Vector Addition Systems and Petri Nets. A k -dimensional *labeled vector addition system*, or *labeled VAS* $M = (A, T, \ell, s, t)$ over alphabet A consists of a set of *transitions* $T \subseteq \mathbb{Z}^k$, a labeling $\ell : T \rightarrow A \cup \{\varepsilon\}$, where ε stands for the empty word and *source* and *target* vectors $s, t \in \mathbb{N}^k$. A labeled VAS defines a transition relation on the set \mathbb{N}^k of *markings*. For two markings $u, v \in \mathbb{N}^k$ we write $u \xrightarrow{a} v$ if there is $r \in T$ such that $u + r = v$ and $\ell(r) = a$, where the addition of vectors is defined as an addition on every coordinate. For two markings $u, v \in \mathbb{N}^k$ we say that u *reaches* v *via a word* w if there is a sequence of markings $u_0 = u, u_1, \dots, u_{n-1}, u_n = v$ such that $u_i \xrightarrow{a_i} u_{i+1}$ for all $i \in \{0, \dots, n-1\}$ and $w = a_0 \cdots a_{n-1}$. For a given labeled VAS M the *language* of M , denoted $L(M)$, is the set of all words $w \in A^*$ such that source reaches target via w . We note that languages of labeled VASs are the same as languages of labeled Petri nets.

Since labeled VAS languages are known to be a full trio (Jantzen, 1979), we only need to prove decidability of the diagonal problem. For instance, this can be done by using the computability of downward closures for VAS languages (Habermehl et al., 2010).

Here, we present an alternative approach to this problem. First we will show that it is enough to consider VASs in which the target marking equals $(0, \dots, 0)$. To this end, let M be a k -dimensional labeled VAS with source vector $s = (s_1, \dots, s_k)$ and target vector $t = (t_1, \dots, t_k)$. We transform M to a new VAS M' in which we add two auxiliary coordinates, called *life* coordinates. The source coordinate is enriched by 0 on one life coordinate and by 1 on the other one, so it is $s' = (s_1, \dots, s_k, 0, 1) \in \mathbb{N}^{k+2}$. Every original transition has two copies. One of these transitions subtracts one from the first life coordinate and adds one to the second life coordinate, the second transition does the opposite. Note that nonemptiness of life coordinates serve just as a necessary condition for firing any transition, as every transition subtracts one from one of these coordinates. Therefore, the original source marking s reaches the original target marking t via the same set of words by which the new source marking s' reaches either $(t_1, \dots, t_k, 0, 1)$ or $(t_1, \dots, t_k, 1, 0)$. We add also two *final* transitions, which subtract the original target vector, subtract one from one of the life coordinates and are labeled ε . Therefore, s can reach t by a word w in M if and only if s' can reach 0^{k+2} by w in M' . Indeed, the implication from left to right is immediate. On the other hand, in order to reach the marking 0^{k+2} in M' , the last transition has to be the final transition, so implication from right to left also holds. Thus it is enough to solve the diagonal problem for VASs in which the target marking is $(0, \dots, 0)$.

We show that this diagonal problem is decidable by a reduction to the place-boundedness problem for VASs with one zero test, which is decidable due to Bonnet et al. (Bonnet et al., 2012). We modify the considered VAS in the following way. For every symbol $a \in A$ we add a new *symbol-coordinate*, which is counting how many times we read symbol a . That is, for every transition which is labeled by $a \in A$, we add 1 in the symbol-coordinate corresponding to a and 0 in the symbol-coordinates corresponding to other symbols. The set of symbol-coordinates computes the Parikh image of a word. We also add one new *minimum-coordinate* and a new transition, labeled by ε , which subtracts one from all the symbol-coordinates and adds one to the minimum-coordinate. It is easy to see that minimum-coordinate can maximally reach the minimum number from the Parikh image tuple. Additionally, for every symbol-coordinate, we add a transition labeled by ε , which can decrease this coordinate by one. The diagonal problem for the original VAS language is equivalent to the question of whether for infinitely many $n \geq 0$ the source marking, enriched by zeros in the new coordinates, reaches a marking $(0, \dots, 0, n)$, with zeros everywhere beside the minimum-coordinate with number n . This can be easily reduced to the place-boundedness for a VAS with one zero test: to check whether there are zeros everywhere else than the minimum-coordinate, one adds yet another coordinate that stores the sum of the coordinates to be checked to 0. Under this condition, it remains to check that the minimum-coordinate can get unbounded. This finishes the proof of decidability of the diagonal problem for languages of labeled VASs.

Effective Parikh equivalence. We mention another obvious consequence of our characterization. Assume separability by PTL is decidable for a full trio \mathcal{D} . Furthermore, suppose that for each given language L in a full trio \mathcal{C} , one can effectively produce a Parikh equivalent language in \mathcal{D} . Then, separability by PTL is also decidable for \mathcal{C} . For example, this means that separability by PTL is decidable for *matrix languages* (Dassow and Păun, 1989), a natural language class that generalizes VAS languages and context-free languages. Indeed, it is well-known that given a matrix language, one can construct a Parikh equivalent VAS language (Dassow and Păun, 1989).

Mixed instances. Our result yields further that if separability by PTL is decidable for each of the full trios \mathcal{C} and \mathcal{D} , then it is also decidable whether given $K \in \mathcal{C}$ and $L \in \mathcal{D}$ are separable by a PTL. Indeed, if the diagonal problem is decidable for \mathcal{C} and for \mathcal{D} , then combining the respective algorithms yields decidability of the diagonal problem for the full trio $\mathcal{C} \cup \mathcal{D}$. For instance, separability of a context-free language from the language of a labeled vector addition system is decidable.

Corollary 5.1. *Let \mathcal{C} and \mathcal{D} be full trios. If separability by PTL is decidable for \mathcal{C} and for \mathcal{D} , then it is also decidable for $\mathcal{C} \cup \mathcal{D}$.*

6. UNDECIDABLE CLASSES

Given the fact that separability by PTL is decidable for context-free languages, the question arises whether the same is true of the context-sensitive languages. Here we show that this is not the case, because of the following reasons. First of all, context-free languages are closed under complementation. Therefore, decidable separability of context-free languages by PTL would imply that it is decidable to test if a given context-sensitive language is PTL. The latter question, however, is already undecidable for context-free languages (Corollary 6.3). In this section we provide these observations with more detail and also exhibit a more general technique to show undecidable separability by PTL.

Thomas Place pointed out to us that, by a proof that follows similar lines the proof of Greibach's Theorem (Greibach, 1968), one can show that testing whether a given context-free language is piecewise testable is undecidable (Place, 2015a). We give the proof below (Proposition 6.2).

We noticed that the proof he gave us even applies to all full trios that contain the language $L_{ab} = \{a^n b^n \mid n \geq 0\}$. Note that this means that for *all the language classes mentioned in Section 5*, it is undecidable whether a given language is piecewise testable.

We introduce some additional material to allow for this generalization. Given a language L , the *full trio generated by L* is the smallest full trio containing L and we denote it by $\hat{\mathcal{M}}(L)$. Since the intersection of full trios is a full trio and every full trio includes the regular languages, such a class indeed exists. Observe that if $L \neq \emptyset$, then $\hat{\mathcal{M}}(L)$ consists of precisely those languages of the form RL , where R is a rational transduction. In particular, a language RL in $\hat{\mathcal{M}}(L)$ can be denoted using a representation for R . Furthermore, note that $\hat{\mathcal{M}}(L)$ is always closed under union, because $RL \cup SL = (R \cup S)L$ and the union of rational transductions is again a rational transduction.

Lemma 6.1. *The universality problem is undecidable for $\hat{\mathcal{M}}(L_{ab})$.*

Proof. We reduce Post's Correspondence Problem (PCP) to the universality problem of $\hat{\mathcal{M}}(L_{ab})$. An instance of PCP consists of an alphabet A and two homomorphisms $\alpha, \beta: A^* \rightarrow \{0, 1\}^*$ and the problem then asks whether there is a $w \in A^+$ with $\alpha(w) = \beta(w)$.

Suppose we are given a PCP instance with $A \cap \{0, 1\} = \emptyset$. Given any homomorphism $\gamma: A^* \rightarrow \{0, 1\}^*$, we first show that one can construct a rational transduction R_γ with

$$R_\gamma L_{ab} = \{uv \mid u \in A^+, v \in \{0, 1\}^*, v \neq \gamma(u)\}.$$

The intuition is that one uses the input word $a^n b^n$ to produce all words uv where the n -th position of v differs from the n -th position of $\gamma(u)$. More precisely, the transducer non-deterministically produces words u and v such that $\gamma(u) \neq v$. Observe that $\gamma(u) \neq v$ means that

- (1) either $|\gamma(u)| > |v|$,
- (2) or $|\gamma(u)| < |v|$,
- (3) or $\gamma(u)$ and v differ at some common position.

Thus, the transducer R_γ is the union of three transducers. We illustrate its construction for the third case. The transducer outputs $u = xyz$ in three steps:

- It outputs some word $x \in A^*$ and, at the same time, reads $a^{|\gamma(x)|}$.
- It outputs a symbol $y \in A$ and remembers in its state a position k in $\gamma(y)$ and the symbol i in $\gamma(y)$ at position k .
- It outputs some word $z \in A^*$.

The transducer produces $v = rst$ in three steps:

- It outputs some word $r \in \{0, 1\}^*$ and, at the same time, reads $b^{|r|}$.
- It outputs a word s that, at position k , has a symbol that differs from i .
- It outputs some word $t \in \{0, 1\}^*$.

Since the input is always $a^n b^n$, the above procedure assures that $|\gamma(x)| = |r|$ and therefore produces precisely the words uv with $v \neq \gamma(u)$. This means that R_γ can be constructed from α and β .

Let $B = A \cup \{0, 1\}$ and $K = B^* \setminus (A^+ \{0, 1\}^*)$. Since $\hat{\mathcal{M}}(L_{ab})$ is effectively closed under union and includes the regular languages, it effectively contains

$$M = R_\alpha L_{ab} \cup R_\beta L_{ab} \cup K.$$

Now clearly, $M = B^*$ if and only if there is no $w \in A^+$ with $\alpha(w) = \beta(w)$. □

Proposition 6.2 (Place, 2015a). *Let \mathcal{C} be a full trio that contains the language L_{ab} . Then, testing whether a given language from \mathcal{C} is piecewise testable is undecidable.*

Proof. Since $\hat{\mathcal{M}}(L_{ab})$ is effectively contained in \mathcal{C} , it clearly suffices to show that piecewise testability is undecidable for $\hat{\mathcal{M}}(L_{ab})$.

Let L be a language from $\hat{\mathcal{M}}(L_{ab})$ over an alphabet A . We show that, if we can decide whether L is piecewise testable, then we can also decide whether L is universal. Since universality is undecidable for $\hat{\mathcal{M}}(L_{ab})$ by Lemma 6.1, we have a contradiction.

Fix K as any language in $\hat{\mathcal{M}}(L_{ab})$ that is not piecewise testable and $\#$ as a symbol that is not in A . We claim that the language $L' = K\#A^* \cup A^*\#L$ is piecewise testable if and only if L is universal. Note that L' belongs to $\hat{\mathcal{M}}(L_{ab})$ since the latter is effectively closed under union.

If L is universal, then $L' = A^*\#A^*$, which is trivially piecewise testable.

If L is not universal, assume by contradiction that L' is piecewise testable. By hypothesis there is a word $w \notin L$. It is then immediate that $K = L'(\#w)^{-1}$ is also piecewise testable (as piecewise testable languages are closed under residuals) which is a contradiction by choice of K . □

Since the context-free languages are a full trio and contain L_{ab} , we have the following as one example. However, note that this is also true for all language classes mentioned in Section 5.

Corollary 6.3. *Testing if a context-free language is piecewise testable is undecidable.*

7. CONCLUDING REMARKS

Since the decidability results we presented seem to be in strong contrast with the remark of Hunt in the introduction, we briefly comment on this. What we essentially do is show that undecidable emptiness-of-intersection for a class \mathcal{C} does not always imply undecidability for separability of \mathcal{C} with respect to some nontrivial class of languages. In the case of separability with respect to piecewise

testable languages, the main reason is basically that we only need to construct intersections of languages from \mathcal{C} with languages that are regular (or even piecewise testable). Here, the fact that such intersections can be effectively constructed, together with decidable emptiness and diagonal problems seem to be sufficient for decidability.

Regarding future work, we see many interesting directions and new questions. Which language classes have a decidable diagonal problem? Which other characterizations are there for decidable separability by PTL? Can Theorem 2.5 be extended to also give complexity guarantees? Can we find similar characterizations for separability by subclasses of PTL, as considered in (Hofman and Martens, 2015)?

Finally, while we now know that separability of context-free languages by PTL is decidable, we do not know what is the complexity. For example, it is not known if the problem is elementary or not.

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